

Approximation Algorithms for 3D Orthogonal Knapsack

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Revised July 10, 2008.

Abstract We study non-overlapping axis-parallel packings of 3D boxes with profits into a dedicated bigger box where rotation is either forbidden or permitted, and we wish to maximize the total profit. Since this optimization problem is NP-hard, we focus on approximation algorithms. We obtain fast and simple algorithms for the non-rotational scenario with approximation ratios $9 + \epsilon$ and $8 + \epsilon$, as well as an algorithm with approximation ratio $7 + \epsilon$ that uses more sophisticated techniques; these are the smallest approximation ratios known for this problem. Furthermore, we show how the used techniques can be adapted to the case where rotation by 90° either around the z -axis or around all axes is permitted, where we obtain algorithms with approximation ratios $6 + \epsilon$ and $5 + \epsilon$, respectively. Finally our methods yield a 3D generalization of a packability criterion and a strip packing algorithm with absolute approximation ratio $29/4$, improving the previously best known result of $45/4$.

Keywords approximation algorithm, computational and structural complexity, geometric configurations

1 Introduction

Given a list $L = \{R_1, \dots, R_n\}$ of boxes with sizes $R_i = (x_i, y_i, z_i)$ and positive profits p_i for each $i \in \{1, \dots, n\}$ and a dedicated bigger box $Q = (a, b, c)$, we study non-overlapping axis-parallel packings of sublists of L into Q which we call *feasible*. For simplicity we call Q a *bin*. We wish to select a sublist which permits a packing and maximizes the profit. This problem will be called the *orthogonal three-dimensional knapsack problem* or *OKP-3*, for short, and we denote the optimal profit by OPT . It is a natural generalization of the knapsack problem (KP) which is known to be NP-hard. This makes an exact algorithm with a polynomial worst-case runtime bound impossible unless $P = NP$ holds. For this reason, we concentrate on approximation algorithms; we refer the reader to [1] for a detailed description of the approach and common notions. W.l.o.g., we assume $a = b = c = 1$ and that each $R_i \in L$ can be packed by otherwise removing infeasible boxes and scaling in $\mathcal{O}(n)$ time. Note that the scaling is only possible in the non-rotational case. In the rotational case the fixed size of the bin is an explicit assumption.

Related Problems. Different geometrically-constrained two- and three-dimensional packing problems were studied, resulting in three main directions.

In *strip packing*, the target area is a strip of infinite height; the objective is to minimize the height of the packing. For the 2D case, Sleator^[2] proved an approximation ratio of $5/2$; Baker *et al.*^[3] obtained an asymptotic approximation ratio of $5/4$. The best known absolute approximation ratio of 2 was obtained independently with different techniques by Schiermeyer^[4] and Steinberg^[5]. Kenyon and Rémila^[6] found an AFP-TAS (asymptotic fully polynomial time approximation scheme) for the problem, and they obtained an additive error of $\mathcal{O}(1/\epsilon^2)$. This additive error was later improved by Jansen and Solis-Oba in [7], where the authors presented an APTAS (asymptotic polynomial time approximation scheme) with additive error 1. For the 3D case, research has focused mainly on the *asymptotic* approximation ratio, where Miyazawa and Wakabayashi^[8] found an algorithm with asymptotic approximation ratio of $27/10$ at most. An asymptotic ratio of $2 + \epsilon$ was obtained by Jansen and Solis-Oba^[9]; this was improved to 1.691 by Bansal *et al.*^[10] The best known *absolute*

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Research supported in part by DFG Project, Entwicklung und Analyse von Approximativen Algorithmen für Gemischte und Verallgemeinerte Packungs- und Überdeckungsprobleme, JA 612/10-1, in part by the German Academic Exchange Service DAAD, in part by project AEOLUS, under EU Contract No. 015964, and in part by a Grant “DAAD Doktorandenstipendium” of the German Academic Exchange Service DAAD. Part of this work was done in duration of visit to the LIG, Grenoble University.

approximation ratio of $45/4$ follows from an asymptotic approximation ratio of $13/4$ from [11]. Li and Cheng^[12] studied the *on-line* version, resulting in a competitive ratio of $29/10$.

In *bin packing*, the objective is to minimize the number of identical bins. For the 1D case, an APTAS was presented by Fernandez de la Vega and Lueker^[13], while Li and Yue^[14] proved the bound of $11/9\text{OPT} + 7/9$ for the popular FFD (first-fit decreasing height) algorithm. The study of FFD was settled recently by Dósa^[15] who proved that the bound $11/9\text{OPT} + 6/9$ is tight. For the 2D case, an asymptotic approximation ratio of 1.691 was obtained by Caprara^[16]. This result was improved to an asymptotic approximation ratio of 1.525 by Bansal *et al.*^[17] Furthermore, Bansal *et al.*^[18] proved that 2D bin packing does not admit an APTAS (asymptotic polynomial time approximation scheme), and therefore there is no FPTAS (fully polynomial time approximation scheme) if $P \neq NP$. They also presented an APTAS for packing d -dimensional cubes into the minimum number of unit cubes in the same paper.

In *knapsack* scenario, the number of bins is a fixed constant^[19], usually 1. For the 2D case, Jansen and Zhang^[20] obtained an approximation ratio of $2 + \epsilon$. Classical 1D knapsack problems are relatively well understood, see [21], [22] for surveys.

Although these problems are closely related, the results cannot be transferred directly. One main contrast between bin and strip packing on the one hand and knapsack on the other hand is that in the first setting all boxes of the instance must be packed but in the latter a feasible selection of items must be found.

Previous Results and Applications. Harren^[23] obtained a ratio of $9/8 + \epsilon$ for the special case of packing cubes, and Chlebík and Chlebíková proved the APX-completeness of the general case^[24].

A *cutting stock* application is cutting blocks with given profits from larger pieces of material to maximize the profit. Another application is the problem of selecting boxes to be transported in a container. Besides these, the problem is motivated from multiprocessor scheduling on grid topology. In this perspective, for a time slice of fixed duration, a set of jobs to be executed must be chosen and each job requires a subgrid of prespecified rectangular shape. For a special case of this application, Ye and Zhang^[25] presented an on-line algorithm, see [26] for a study of similar problems.

New Results. Our contribution is a fast and simple $(9 + \epsilon)$ -approximation algorithm based on strip packing which is refined to an $(8 + \epsilon)$ -approximation algorithm. Both of these have practical running times. With more sophisticated techniques, we obtain a $(7 + \epsilon)$ -approximation

algorithm. We also study the case where rotation by 90° either around the z -axis or around all axes is permitted, where we significantly improve upon the approximation ratios of the algorithms presented for OKP-3. We derived approximation ratios of $6 + \epsilon$ for the former and $5 + \epsilon$ for the latter case. Finally our methods yield a three-dimensional generalization of a packability criterion and a strip packing algorithm with absolute approximation ratio $29/4$.

This paper is organized as follows. In Section 2, we present a fast algorithm for non-rotational packing which is improved in Section 3. In Section 4 we generalize a two-dimensional packability criterion for boxes, and improve a known result on strip packing in Section 5 before turning back to the knapsack problem in Section 6, where we obtain a better yet more costly algorithm. Finally, we discuss the cases of rotational packing in Section 7, and conclude the paper with open problems in Section 8.

2 Algorithm Based on Strip Packing

We approximately solve a relaxation by selecting $L' \subseteq L$ that is at least near-optimal and has a total volume of at most 1. This relaxed solution is partitioned into 9 sublists. For each of these a packing into the bin will be generated. Out of these, one with the maximum profit is chosen, resulting in a $(9 + \epsilon)$ -approximation algorithm. More precisely L' will be packed into a strip $[0, 1] \times [0, 1] \times [0, \infty)$ by a level-oriented algorithm, i.e., an algorithm which packs all boxes into disjoint levels and stacks these levels on top of one another into the strip. We partition the strip into packings of sublists of L' and among these return one with the maximum profit.

For each box R_i , the rectangle (x_i, y_i) is called the *base rectangle* of R_i , denoted as $br(R_i)$. Such a rectangle (x_i, y_i) is called

$$\begin{aligned} \textit{big} &: \Leftrightarrow x_i \in (1/2, 1] \quad \text{and} \quad y_i \in (1/2, 1], \\ \textit{long} &: \Leftrightarrow x_i \in (1/2, 1] \quad \text{and} \quad y_i \in (0, 1/2], \\ \textit{wide} &: \Leftrightarrow x_i \in (0, 1/2] \quad \text{and} \quad y_i \in (1/2, 1], \\ \textit{small} &: \Leftrightarrow x_i \in (0, 1/2] \quad \text{and} \quad y_i \in (0, 1/2]. \end{aligned}$$

For each list L of boxes we use $V(L)$ to denote the total volume of L , and for each list \hat{L} of rectangles use $A(\hat{L})$ to denote the total area of \hat{L} . Furthermore, $P(L)$ denotes the total profit of L . Finally, for each list L of boxes we use $H(L)$ to denote the height of a packing of L where the packing itself will be clear from the context. We use the following theorem from [20] which is a refinement of the main result from [5].

Theorem 1^[20]. *Let L be a list of n rectangles such that $A(L) \leq 1/2$ holds, and either no long or no wide rectangles occur in L , i.e., there might be either long or wide rectangles, but not both. Then L permits a feasible packing into the unit square which can be generated in time $\mathcal{O}(n \log^2 n / \log \log n)$.*

Note that rectangles which are long or wide are not big. It is therefore possible in Theorem 1 that there is a big rectangle in L . First we apply the modified strip packing algorithm, and then we construct the partition of the strip. The strip packing algorithm uses Theorem 1 to obtain an *area guarantee* for each level but the last, improving a result from [11].

Algorithm A

1. Partition L into $L_1 := \{R_i \mid \text{br}(R_i) \text{ is long}\}$ and $L_2 := L \setminus L_1$. W.l.o.g., we have

$$L_1 = \{R_1, \dots, R_m\}$$

and

$$L_2 = \{R_{m+1}, \dots, R_n\}.$$

2. Generate the packing for L_1 as follows.

- 2.1. Find the boxes R_i in L_1 for which the area of $\text{br}(R_i)$ is greater than $1/4$, which are R_{p+1}, \dots, R_m , w.l.o.g., Stack these on top of one another in direction z , each on its own level.
- 2.2. Sort the remaining boxes R_1, \dots, R_p in non-increasing order of z_i , resulting in a list L'_1 .
- 2.3. Partition L'_1 into consecutive sublists L''_1, \dots, L''_v where the total base area of each sublist is as close to $1/2$ as possible but not greater. Pack each of these sublists on a level by itself using Theorem 1. Stack all of these levels on top of one another in direction z .

3. Generate the packing for L_2 in a similar way as for L_1 by Theorem 1. The resulting steps are called Steps 3.1~3.3.

4. Concatenate the packings of L_1 and L_2 to obtain a packing of L .

Theorem 2. *For each list L of n boxes, Algorithm A generates a packing of height at most $4V(L) + Z_1 + Z_2$, where Z_1 and Z_2 are the heights of the first levels generated in Steps 2.3 and 3.3. The construction can be carried out in time $\mathcal{O}(n \log^2 n / \log \log n)$.*

Proof. Consider Step 2.3 which generates levels L''_1, \dots, L''_v . Let h_i denote the height of level L''_i for each $i \in \{1, \dots, v\}$. Each box R_i for $i \in \{p+1, \dots, m\}$ processed in Step 2.1 forms a level itself, so the total height for the packing of L_1 is

$$H(L_1) := \sum_{i=p+1}^m z_i + \sum_{i=1}^v h_i, \quad (1)$$

and we have $V(L) = V(L_1) + V(L_2)$. Furthermore,

$$V(L_1) = \sum_{i=p+1}^m x_i y_i z_i + \sum_{i=1}^v V(L''_i)$$

holds. In Step 2.1 we assert that $x_i y_i > 1/4$ for $i \in \{p+1, \dots, m\}$, so we get

$$\sum_{i=p+1}^m x_i y_i z_i > \frac{1}{4} \sum_{i=p+1}^m z_i. \quad (2)$$

For each $i \in \{1, \dots, v-1\}$, the total base area of the boxes in L''_i is larger than $1/4$ since otherwise the next box would have also been put in L''_i in Step 2.3. As the boxes are sorted in non-increasing order of height, and thus $z_j \geq h_{i+1}$ for $R_j \in L''_i$, we get that

$$\sum_{R_j \in L''_i} x_j y_j > \frac{1}{4} \quad (3)$$

holds, hence

$$V(L''_i) \geq \frac{1}{4} h_{i+1} \quad (4)$$

for each $i \in \{1, \dots, v-1\}$. Combining (2) and (4) results in

$$V(L_1) > \frac{1}{4} \sum_{i=p+1}^m z_i + \frac{1}{4} \sum_{i=2}^v h_i$$

which together with (1) implies that

$$H(L_1) < 4V(L_1) + h_1 \leq 4V(L_1) + Z_1 \quad (5)$$

holds. For Step 3 we get

$$H(L_2) < 4V(L_2) + Z_2 \quad (6)$$

with a very similar analysis. From (5) and (6) we conclude that the height of our packing of L is bounded from above by

$$H(L_1) + H(L_2) < 4V(L) + Z_1 + Z_2.$$

The runtime is dominated by the application of Theorem 1 and thus bounded by $\mathcal{O}(n \log^2 n / \log \log n)$. \square

The second part of the algorithm is a partition of the arrangement generated by Algorithm A into at most 9 bins, see Fig.1.

Algorithm B

1. Set $\delta := \epsilon/(9 + \epsilon)$. Use an FPTAS for KP from [21, 27] to select $L' \subseteq L$ such that $V(L') \leq 1$ and $P(L') \geq (1 - \delta)OPT$ holds, where OPT denotes the optimum of the generated KP instance.
2. Use Algorithm A to generate a packing of L' into the strip but separate the first levels generated in Steps 2.3 and 3.3. Pack these into a bin each.

3. By Theorem 2 the remaining strip has a height of at most $4V(L') \leq 4$. Consider the cutting unit squares $[0, 1] \times [0, 1] \times \{i\}$ for $i \in \{1, 2, 3\}$. Generate a partition of the region $[0, 1] \times [0, 1] \times [0, 4]$ into 7 subsets, i.e., 4 subsets which are positioned each in the regions $[0, 1] \times [0, 1] \times [i-1, i]$ for $i \in \{1, \dots, 4\}$ but not intersecting any of the unit squares, and 3 subsets of boxes each of which intersects with one of the three cutting unit squares.
4. Out of the sets generated in Steps 2 and 3 returns one with the maximum profit.

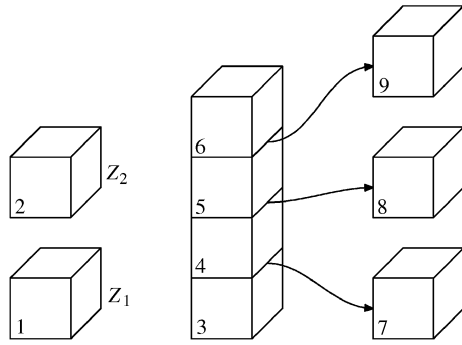


Fig.1. At most 9 bins are generated by Algorithm B.

Each set generated in Steps 2 and 3 permits a feasible packing into the unit cube which is available as a byproduct of Algorithm A. L' is partitioned into at most 9 subsets by Algorithm B, as illustrated in Fig.1.

Theorem 3. *Algorithm B is a $(9 + \epsilon)$ -approximation algorithm for OKP-3 with running time $\mathcal{O}(T_{KP}(n, \epsilon) + n \log^2 n / \log \log n)$, where $T_{KP}(n, \epsilon)$ is the running time of the FPTAS used for KP from [21, 27]; furthermore, this bound is tight. Note that in [21] the best running time for an FPTAS for the knapsack problem is*

$$\mathcal{O}(n \min\{\log n, \log(1/\epsilon)\} + 1/\epsilon^2 \log(1/\epsilon) \min\{n, 1/\epsilon \log(1/\epsilon)\}).$$

Proof. Clearly $9 + \epsilon$ is an upper bound for the ratio and the running time is dominated by solving the knapsack instance and by Algorithm A since

$$\frac{(1 - \delta)}{9} = \frac{1}{9 + \epsilon}$$

holds. For the following instance this bound can be attained. We have 10 boxes $R_1 := (1/2, 1/2, 2/15)$, $R_2 := (1, 1/4, 2/15)$, $R_3 := (1, 2/7, 3/4)$, $R_4 := \dots := R_7 := (1, 2/7, 1/2)$, $R_8 := (1, 2/7, 1/4 + 2/15)$, $R_9 := (1, 2/7, 2/15)$ and $R_{10} := (1, 1, 1)$. Furthermore, $p_1 := \dots := p_9 := 1/(9 + \epsilon)$, and $p_{10} := 1$. Let $S_1 := \{R_1, \dots, R_9\}$ and $S_2 := \{R_{10}\}$. It is clear that

S_2 is an optimal solution. Elementary calculation shows $V(S_1) = 1$ and $P(S_1) = 1 - \delta$, and hence S_1 may be selected in Step 1 of Algorithm B. Application of Algorithm B, assuming that the boxes are stacked in increasing order of index in Step 2.1 of Algorithm A, yields 9 bins each containing an item with profit $1/(9 + \epsilon)$, see Fig.2. \square

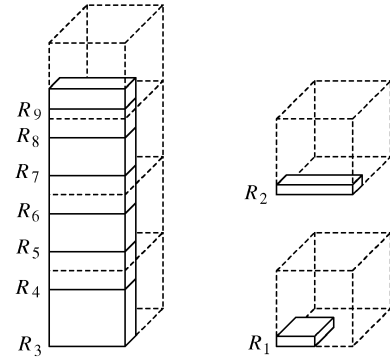


Fig.2. Packing of S_1 with Algorithm B.

Note that only the subset that is returned needs to be packed level-wise using the algorithm from Theorem 1 while the discarded subsets need not be arranged. Algorithm B can be used to solve the special cases where we wish to maximize the number of selected boxes or the volume by setting $p_i := 1$ or $p_i := x_i y_i z_i$ for each $i \in \{1, \dots, n\}$. This also holds for the other algorithms which we present.

In [21, 27] approximation algorithms for various knapsack problems are found. Using them, Algorithm B can be generalized by replacing the KP solver in Step 1, yielding algorithms for *unbounded* OKP-3 and *multiple-choice* OKP-3, see [21, 27] for notions and details. Algorithm B can be modified to yield a ratio of 18 with a much better running time by using a 2-approximation algorithm for classical KP from [21, 27], thus replacing $T_{KP}(n, \epsilon)$ by $\mathcal{O}(n)$ in Theorem 3.

3 Refined Construction

In this section we show how to refine Algorithm B to yield an approximation ratio of $(8 + \epsilon)$. We identify two possible improvements on Algorithm A: first, by increasing the area guarantee and second, by decreasing the heights Z_1 and Z_2 of the additional strip packing levels.

In Algorithm A and the proof of Theorem 2, the area bound $1/2$ from Theorem 1 was used. We separated boxes with base area greater than $1/4$, resulting in the area guarantee of $1/2 - 1/4 = 1/4$ for each level generated in Steps 2.2 and 2.3 except the last one. By

improving the area guarantee we will improve the height bound of the strip.

So far we arbitrarily chose direction z to be the axis for level generation, but any direction $d \in \{x, y, z\}$ will do. This has the advantage that packing in a direction where all boxes are short, i.e., at most $1/2$, implies that the heights of the additional levels are bound by $1/2$.

Let us introduce the notion of big boxes in certain directions: for any direction $d \in \{x, y, z\}$ a box R_i is called d -big: $\Leftrightarrow d_i \in (1/2, 1]$, and we use X, Y and Z to denote the set of boxes that are d -big for the corresponding direction. Any box that is d -big for every direction $d \in \{x, y, z\}$ will be called a *big* box. Finally, a box R_i is called *small*: $\Leftrightarrow x_i \in (0, 1/2], y_i \in (0, 1/2]$, and $z_i \in (0, 1/2]$.

The remainder of the section is organized as follows. First, we give a refined version of Algorithm A that is restricted to packing small boxes. Second, we show how to partition the boxes which are not small into three sets according to the big direction. Third, we give the overall algorithm which is based on the partition into d -big and small boxes.

The following algorithm is applied only on *small* items.

Algorithm C

1. Find the boxes R_i in L for which the area of $br(R_i)$ is greater than $1/10$, and they are R_1, \dots, R_m , w.l.o.g., Sort these in non-increasing order of z_i , resulting in a list L_1 . Arrange these in groups of 4 boxes each, except for the last group. Each group can be put on a separate level by placing the boxes into the corners of the level. Stack these levels on top of one another in direction z .
2. Sort the remaining boxes R_{m+1}, \dots, R_n in non-increasing order of z_i , resulting in a list L_2 .
3. Partition L_2 into consecutive sublists L'_1, \dots, L'_v where the total base area of each sublist is as close to $1/2$ as possible but not greater. Pack each of these sublists on a level by itself using Theorem 1. Stack all of these levels on top of one another in direction z .
4. Concatenate the packings of L_1 and L_2 to obtain a packing of L .

Note that we formed two groups and obtained an area guarantee of $2/5 = 1/2 - 1/10$ for each layer except the last ones generated in Steps 1 and 3. To avoid confusion we point out that the area guarantee does not hold for the *last* generated layers, while the summands Z_1 and Z_2 in Theorem 2 are the heights of the respective *first* layers. Similarly to the proof of Theorem 2, we obtained the following results using the area guarantee of $2/5$.

Theorem 4. *For each list L of n small boxes, Algorithm C generates a feasible packing of height at*

most

$$5/2V(L) + Z_1 + Z_2,$$

where $Z_1 \leq 1/2$ and $Z_2 \leq 1/2$ are the heights of the first levels generated in Steps 1 and 3. The construction can be carried out in time $\mathcal{O}(n \log^2 n / \log \log n)$.

Lemma 5. *Each list L of n small boxes with $V(L) \leq 1$ permits a feasible packing into at most 5 bins. The construction can be carried out in time $\mathcal{O}(n \log^2 n / \log \log n)$; the bound of 5 is tight for the used construction.*

Proof. Use Algorithm C to arrange L in a strip, but separate the first levels generated in Steps 1 and 3. Since L contains only small boxes, these two levels can be packed together into a bin. By Theorem 4, the remaining strip has a height of at most $5/2$. Consider the two cutting unit squares $[0, 1] \times [0, 1] \times \{i\}$ for $i \in \{1, 2\}$. Generate a partition of the region $[0, 1] \times [0, 1] \times [0, 5/2]$ into 5 subsets, namely first 3 subsets which are each positioned in the regions $[0, 1] \times [0, 1] \times [i-1, i]$ for $i \in \{1, 2\}$ as well as the region $[0, 1] \times [0, 1] \times [2, 5/2]$ but not intersecting any of the unit squares, and furthermore the rest 2 subsets of boxes each of which intersects with one of the two cutting unit squares. The first three sets can be packed into one bin each. Since L contains only small boxes, the last two sets can be arranged together into one additional bin by aligning them at the top and bottom of the bin, respectively. We have at most 5 bins, see Fig.3. The running time is dominated by Algorithm C and thus bounded by $\mathcal{O}(n \log^2 n / \log \log n)$.

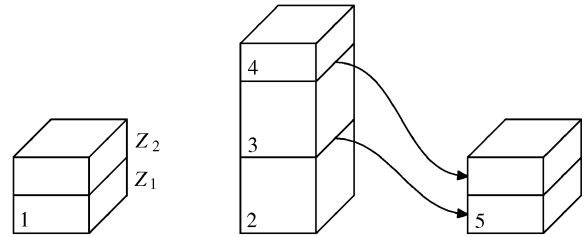


Fig.3. Small boxes can be packed into at most 5 bins.

To show the tightness of the bound, let $\gamma := 1/500$ and consider the instance L consisting of $R_1 := \dots := R_{29} := (1/2, 1/5 + \gamma, 1/3 + \gamma)$, $R_{30} := (\gamma, \gamma, 1/2)$, $R_{31} := \dots := R_{33} := (1/2, 1/5, \gamma)$ and $R_{34} := (1/2, 1/5 - 2\gamma^2, \gamma)$. Note that

$$\begin{aligned} V(L) &= 29/30 + 122/15\gamma + 15\gamma^2 - \gamma^3 \\ &< 29/30 + 1/60. \end{aligned}$$

Application of Algorithm C packs R_1, \dots, R_{29} in Step 1, resulting in 8 layers. All remaining boxes are packed into one more layer in Step 3. The height of each layer

is greater than $1/3$, which means that the layers cannot be arranged in less than 5 bins. \square

A partition of the boxes which are not small is given by the following lemma. In the sequel we use the notion of the *projection* of a box; for each box $R = (x, y, z)$ we call the rectangle (y, z) the x -projection, the rectangle (x, z) the y -projection, and the rectangle (x, y) the z -projection of R .

Lemma 6. *Let L be a list of n boxes in which no small boxes and at most 3 big boxes occur. Then L can be partitioned into sets X' , Y' and Z' in time $\mathcal{O}(n)$, such that each of these contains at most one big box, and the x -projections of boxes in X' , the y -projections of boxes in Y' and the z -projections of boxes in Z' contain no long or no wide rectangles.*

Proof. Remove the at most 3 big boxes from L and distribute them in X' , Y' and Z' such that in each of these sets at most one big box occurs. Set $X' := X' \cup \{R_i \in L \mid x_i > 1/2, z_i \leq 1/2\}$, $Y' := Y' \cup \{R_i \in L \mid y_i > 1/2, x_i \leq 1/2\}$ and finally $Z' := Z' \cup \{R_i \in L \mid z_i > 1/2, y_i \leq 1/2\}$ to obtain the claim. To see that X' , Y' and Z' form a partition, consider w.l.o.g., a box R that is x -big. If this box is y -big and z -big, it is big and therefore included in one of the sets. Otherwise it is either z -big and in Z' or small in direction z and thus in X' . On the other hand, no box can be in more than one of the sets. \square

We are now ready to give the overall algorithm. To avoid repetition, we enumerate the cases in the analysis only.

Algorithm D

1. Set $\delta := \epsilon/(8 + \epsilon)$. Use a PTAS for non-geometric 4D KP from [21, 27] to select $L' \subseteq L$ such that $P(L') \geq (1 - \delta)OPT$, where OPT denotes the optimum of the integral linear program

$$\text{maximize } \sum_{i=1}^n p_i R_i \quad \text{subject to } R \in P,$$

where R_i is an indicator variable for the box of the same name. Furthermore we define the set P by the constraints

$$\begin{aligned} \sum_{i=1}^n x_i y_i z_i R_i &\leq 1, \\ \sum_{R_i \in X} y_i z_i R_i &\leq 1, \\ \sum_{R_i \in Y} x_i z_i R_i &\leq 1, \\ \sum_{R_i \in Z} x_i y_i R_i &\leq 1. \end{aligned}$$

In total, P is a polytope of nonnegative integers.

2. Partition L' into at most 8 subsets which permit a feasible packing as described below. Out of these, returns one with the maximum profit.

Theorem 7. *Algorithm D is a $(8 + \epsilon)$ -approximation algorithm for OKP-3 with running time $\mathcal{O}(T_{4DKP}(n, \epsilon) + n \log^2 n / \log \log n)$, where $T_{4DKP}(n, \epsilon)$ is the running time of the PTAS used for 4D KP from [21, 27]; furthermore this bound is tight. Note that in [21] the best running time for a PTAS for the multidimensional knapsack problem is $\mathcal{O}(n^{\lceil d/\epsilon \rceil - d})$.*

Proof. The first constraint of the integral linear program models the volume bound of the box. The other constraints are area bounds for d -big boxes for $d \in \{x, y, z\}$, motivated by the observation that the d -projections of d -big boxes do not overlap. Thus the given program is a relaxation of our problem.

We have not imposed a bound on the number of big boxes in the relaxation, but due to the area conditions there are at most 3 big boxes in the selected set. We consider two cases according to the total projection area of the d -big boxes.

Case 1. There is a direction $d \in \{x, y, z\}$ such that the total d -projection area of all d -big boxes in L' is larger than or equal to $1/2$. In this case all d -big boxes can be packed into at most 3 bins with a construction from [20], which can be done in time $\mathcal{O}(n \log^2 n / \log \log n)$, resulting in a volume of at least $1/4$ being packed. The total volume of the remaining boxes is bounded by $3/4$ and each remaining box has a d -height of at most $1/2$. We apply Algorithm A in direction d which results in a strip of d -height at most 3 and two additional levels of d -height at most $1/2$ each. Similar to the proof of Lemma 5 all these sets can be packed into at most 5 bins, see Fig.3, generating at most 8 bins in total.

Case 2. For each $d \in \{x, y, z\}$ the total projection area of all d -big boxes is smaller than $1/2$. By Lemma 6 we partition the set $\{R_i \in L' \mid R_i \text{ is not small}\}$ into sets X' , Y' and Z' such that the total projection area of X' , Y' and Z' for the corresponding direction is not greater than $1/2$, and the x -projections of boxes in X' , the y -projections of boxes in Y' and the z -projection of boxes in Z' contain no long or no wide rectangles, respectively. Furthermore, each of these sets contains at most one big box. By Theorem 1 the sets X' , Y' and Z' can be packed into at most one bin each, resulting in at most 3 bins in total. Let S denote the set of small boxes; these are not yet packed. Clearly $V(S) \leq 1$ holds, so by Lemma 5 the set S can be packed into at most 5 bins, which results in at most 8 bins in total. The runtime bound follows from the fact that we can distinguish between the two cases in time $\mathcal{O}(n)$.

For the tightness of the bound, consider the instance L in which R_1, \dots, R_{34} are as in the proof of Lemma 5, $R_{35} := (1, 1, 1/180)$, $R_{36} := (1, 1/180, 1)$, $R_{37} := (1/180, 1, 1)$, and $R_{38} := (1, 1, 1)$. The profits are defined by $p_i := 1/[9(8 + \epsilon)]$ for $i \in \{1, \dots, 4, 30, \dots, 34\}$, $p_i := 1/[8(8 + \epsilon)]$ for $i \in \{5, \dots, 28\}$, $p_i := 1/(8 + \epsilon)$ for $i \in \{29, 35, 36, 37\}$, and $p_{38} := 1$. Let $S_1 := L \setminus \{R_{38}\}$ and $S_2 := \{R_{38}\}$. Since $P(S_1) = 8/(8 + \epsilon) = (1 - \delta) < 1 = P(S_2)$, S_2 is an optimal solution. Elementary calculation verifies that S_1 may be chosen in Step 1 of Algorithm D. Application of Algorithm D leads to Case 2 in the analysis above, where $X' = \{R_{35}\}$, $Y' = \{R_{37}\}$ and $Z' = \{R_{36}\}$. Each of these sets is packed into a separate bin. The remaining boxes are small and are packed into 5 bins as in the proof of Lemma 5. In total, 8 bins are generated; the profits are chosen such that each bin yields a profit of exactly $1/(8 + \epsilon)$. \square

4 Packability Criterion

Theorem 1 has a number of applications besides the use in this paper, see [20, 28]. In this section we give a generalization to Theorem 1 for the three-dimensional case.

Lemma 8. *Let L be a list of n boxes such that $V(L) \leq 1/8$ holds, and no d_1 -big and no d_2 -big boxes occur in L for $d_1, d_2 \in \{x, y, z\}$ and $d_1 \neq d_2$. Then L permits a feasible packing into the unit bin which can be generated in time $\mathcal{O}(n \log^2 n / \log \log n)$.*

Observe that similarly to Theorem 1 we are allowed to have items that are d -big in a certain direction d but not in any other direction. A difference between the original two-dimensional version and our generalization is that in the original version a single big item—which is neither long nor wide—was allowed, but is forbidden here.

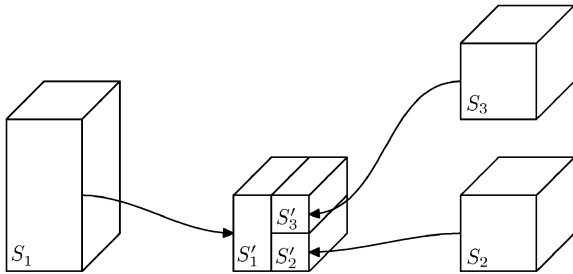


Fig.4. Arrangement of the stacks into a cube.

Proof. Assume w.l.o.g., assume that there are no x -big and no z -big boxes in L . Obtain the set L_{scaled} by scaling the given instance by 2 in directions x and z . We use Algorithm A in direction z to pack the

boxes in L_{scaled} into one stack S_1 of height at most $4V(L_{\text{scaled}}) \leq 4 \cdot 1/2 = 2$ and two additional stacks S_2 and S_3 of height limited by 1 which correspond to the summands Z_1 and Z_2 in Theorem 4. Rescaling everything yields a stack S'_1 of height at most 1 and two additional stacks S'_2 and S'_3 of heights $1/2$. All stacks have a width of at most $1/2$ in direction x . Thus we can arrange all stacks in a unit cube as shown in Fig.4. \square

5 Improved Strip Packing Algorithm

In Theorem 7 in [11], an approximation algorithm for 3D strip packing with asymptotic approximation ratio $13/4$ is presented, and more precisely the bound is $13/4 \text{OPT} + 8Z$, where Z is the height of the highest item. In this section we show that the additive constant of this bound can be easily improved upon by using Theorem 1; more precisely we obtain the following result by using a more suitable subdivision which results in fewer groups.

Theorem 9. *For each list L of n boxes a packing into the strip $[0, 1] \times [0, 1] \times [0, \infty)$ of height at most $13/4 \text{OPT}(L) + 4Z$ can be generated, where OPT denotes the minimum attainable packing height and Z is the height of the highest box. The running time is polynomial in n .*

As we already mentioned in the introduction, the currently best known *asymptotic* approximation ratio for this problem is 1.691 by Bansal et al.^[10] Our result improves the best known *absolute* approximation ratio of $45/4$ which follows from [11] to $29/4$.

Proof. The construction and the proof are similar as in [11] and included for completeness. We partition L into five groups by letting

$$\begin{aligned} L_1 &:= \{R_i \mid \text{br}(R_i) \text{ is big}\}, \\ L_2 &:= \{R_i \mid \text{br}(R_i) \text{ is long and } A(\text{br}(R_i)) \leq 1/6\}, \\ L_3 &:= \{R_i \mid \text{br}(R_i) \text{ is long and } A(\text{br}(R_i)) > 1/6\}, \\ L_4 &:= \{R_i \mid \text{br}(R_i) \text{ is wide or small,} \\ &\quad \text{and } A(\text{br}(R_i)) \leq 1/6\}, \text{ and} \\ L_5 &:= \{R_i \mid \text{br}(R_i) \text{ is wide or small,} \\ &\quad \text{and } A(\text{br}(R_i)) > 1/6\}, \end{aligned}$$

and discuss how to obtain good corresponding area guarantees. Clearly, an area guarantee of $1/4$ can be obtained for L_1 by simply stacking the boxes on top of one another. Furthermore, the groups L_2, \dots, L_5 can be sorted in non-increasing order of z_i . Now for each of these, we proceed similarly as in Step 2.3 of Algorithm A. For L_2 we generate layers using Theorem 1

and obtain an area guarantee of $1/2 - 1/6 = 1/3$ for each layer except the last one. In an even simpler way, the group L_3 can be packed by putting at least two boxes on each layer except for the last one; we obtain an area guarantee of $1/3$ for each but the last layer. By using Theorem 1, we generate layers for the items in L_4 and again obtain an area guarantee of $1/3$ for each layer but the last. Finally, the boxes of L_5 can be packed by placing two of them on each layer except for the last one; again we obtain an area guarantee of $1/3$ for each but the last layer.

Now let $H(L_1)$ denote the height of the packing generated for L_1 and, for each $i \in \{2, \dots, 5\}$, let $H(L_i)$ denote the height of the packing of L_i minus the height of the corresponding *first* layer generated. Let $H(L)$ denote the total height of the packing. For

$$H := \sum_{i=2}^5 H(L_i)$$

clearly

$$H(L) \leq H(L_1) + H + 4Z \quad (7)$$

holds. Furthermore, we have

$$V(L_1) \geq \frac{H(L_1)}{4}.$$

Now let $i \in \{2, \dots, 5\}$ and let L'_1, \dots, L'_j denote the levels generated for L_i ; finally let $V(L'_k)$ denote the total volume of boxes in L'_k and let h_k denote the height of L'_k for each $k \in \{1, \dots, j\}$. Since for L'_1, \dots, L'_{j-1} we have an area guarantee of at least $1/3$, we obtain

$$V(L_i) > \sum_{k=1}^{j-1} V(L'_k) \geq \frac{1}{3} \sum_{k=2}^j h_k = \frac{H(L_i)}{3}.$$

In total, we obtain

$$OPT(L) \geq V(L) = \sum_{i=1}^5 V(L_i) > \frac{H(L_1)}{4} + \frac{H}{3}.$$

Obviously we also have $OPT(L) \geq H(L_1)$, and thus

$$OPT(L) \geq \max \left\{ H(L_1), \frac{H(L_1)}{4} + \frac{H}{3} \right\} \quad (8)$$

holds. Now we study the ratio

$$r := \frac{H(L_1) + H}{\max \left\{ H(L_1), \frac{H(L_1)}{4} + \frac{H}{3} \right\}}. \quad (9)$$

If $H(L_1) \geq H(L_1)/4 + H/3$, which means that $H(L_1) \geq 4/9H$, then we have

$$r = \frac{H(L_1) + H}{H(L_1)} = 1 + \frac{H}{H(L_1)} \leq 1 + \frac{9}{4} = \frac{13}{4}.$$

If $H(L_1) \leq H(L_1)/4 + H/3$, which means that $H(L_1) \leq 4/9H$, then

$$r = \frac{H(L_1) + H}{\frac{H(L_1)}{4} + \frac{H}{3}}$$

holds. In this case, r is a strictly monotonically increasing function of $H(L_1)$, thus the maximum is attained for $H(L_1) = 4/9H$, and $13/4$ is the corresponding maximum value. In total, we have $r \leq 13/4$. Rearrangement of the inequalities above yields

$$\begin{aligned} H(L) &\stackrel{(7)}{\leq} H(L_1) + H + 4Z \\ &\stackrel{(9)}{=} r \max \left\{ H(L_1), \frac{H(L_1)}{4} + \frac{H}{3} \right\} + 4Z \\ &\stackrel{(8)}{\leq} \frac{13}{4} OPT(L) + 4Z \end{aligned}$$

which proves the claim. \square

Since $OPT \geq Z$ we easily derive the following corollary.

Corollary 10. *The absolute approximation ratio of the given algorithm is at most $29/4$.*

6 Enumerations and Shifting Technique

Algorithms B and D generate cutting areas in the strip, resulting in subsets that have to be repacked. We permit further loss of profit by removing more boxes to discard inconvenient layers; the loss will be suitably bounded. The improvement will be at the cost of a considerably larger running time due to a large enumeration; we thus omit a run time analysis. First we introduce a shifting technique to remove sets intersecting the cutting areas and the additional layers. In the sequel we use the notion of a *gap* which is a rectangular region in the strip that does not intersect the interiors of packed boxes.

Lemma 11. *Let $L = \{R_1, \dots, R_n\}$ be a list of boxes with $z_i \leq \epsilon$ for each $R_i \in L$. Suppose L admits a packing into a strip of height at most h and let m be a positive integer. Then we can create m gaps of shape $[0, 1] \times [0, 1] \times [0, \epsilon]$ in a packing of height h by deleting boxes such that for the remaining list $L' \subseteq L$ the inequality $P(L') \geq (1 - 2(m+1)\epsilon/h)P(L)$ holds. The construction can be done in time polynomial in n .*

Proof. Consider the original packing in a strip of height exactly h . We partition the strip into regions

of height ϵ and eventually one region of lower height. More precisely we define $p := \lceil h/\epsilon \rceil$ and partition the strip of height h into p regions S_1, \dots, S_p of shape

$$[0, 1] \times [0, 1] \times [0, \epsilon],$$

where the uppermost region is possibly of lower height. Then for each $i \in \{1, \dots, p\}$ let $T_i = \{R_j \in L \mid R_j \cap S_i \neq \emptyset\}$ and let U_1, \dots, U_{m+1} be the $m+1$ sets out of T_1, \dots, T_p which have the lowest profit. It is easy to see that removing these from the packing causes a loss of profit which is at most $2(m+1)/pP(L)$ since each box is included in at most two of the sets T_1, \dots, T_p . We remove $m+1$ sets, since the uppermost region can be among the sets with the lowest profit, but this region might have height lower than ϵ ; by removing $m+1$ sets we assert that we create at least m gaps of height ϵ . Let L' be the set of remaining boxes; then

$$\begin{aligned} P(L') &\geq P(L) - \frac{2(m+1)}{p}P(L) = \left(1 - \frac{2(m+1)}{p}\right)P(L) \\ &\geq \left(1 - 2(m+1)\frac{\epsilon}{h}\right)P(L) \end{aligned}$$

holds. \square

Note that the construction above can be carried out in any directions.

Theorem 12. *Let R_1, \dots, R_n be a list of boxes with $z_i \leq \epsilon$ for each $R_i \in L$ and $V(L) \leq \alpha \leq 1$. Then it is possible to select $L'' \subseteq L$ such that $P(L'') \geq (1 - 12\epsilon)P(L)$ holds, and L'' admits a feasible packing into at most $\lceil 4\alpha \rceil$ bins. The construction can be carried out in time polynomial in n .*

Proof. W.l.o.g., $\epsilon \leq 1/12$ holds, since otherwise $L'' = \emptyset$ will do. Use Algorithm A to pack L into a strip which is of height $h \leq 4\alpha$ and two additional layers L_1 and L_2 by Theorem 2. If $h \leq 1 - 2\epsilon$ we can clearly pack L into 1 bin without losing any profit. For $1 - 2\epsilon < h < 1$, use Lemma 11 to generate 2 gaps in the strip, which causes a loss of profit of at most $6\epsilon/hP(L) \leq 12\epsilon P(L)$, since $\epsilon \leq 1/12$. Moreover, placing L_1 and L_2 into the gaps gives a feasible packing into 1 bin. For the remainder of the proof we thus assume $h \geq 1$. The following construction is illustrated in Fig.5. Use Lemma 11 to generate at most 5 suitable gaps in the strip, resulting in a loss of profit of at most $12\epsilon/hP(L)$; since $h \geq 1$, this loss is bounded by $12\epsilon P(L)$. The remaining set of boxes in the strip and L_1 and L_2 is denoted as L'' . Consider the 3 cutting unit squares $[0, 1] \times [0, 1] \times \{i\}$ for $i \in \{1, 2, 3\}$, and let L_3, L_4 and L_5 be the sets of boxes in the strip that intersect with these unit squares, respectively. W.l.o.g., none of the sets L_1, \dots, L_5 is empty; otherwise it is removed from consideration. Note that each of the sets

L_1, \dots, L_5 can be arranged on a layer of height at most ϵ , so we generate a feasible packing by arranging them into the 5 gaps. In the resulting packing, the 3 cutting unit squares $[0, 1] \times [0, 1] \times \{i\}$ for $i \in \{1, 2, 3\}$ do not intersect with any box. Furthermore, all layers L_1, \dots, L_5 are merged in the strip; the packing can be rearranged into $\lceil 4\alpha \rceil$ bins. \square

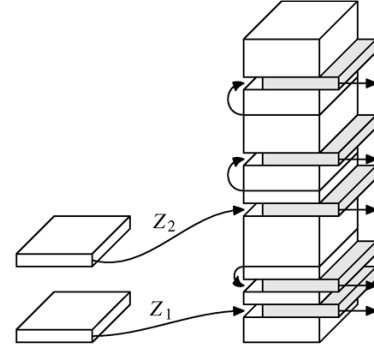


Fig.5. Shifting technique described in Theorem 12.

Like before for any $d \in \{x, y, z\}$ we call a box R_i d - ϵ -big $\Leftrightarrow d_i \in (\epsilon, 1]$ and d - ϵ -small $\Leftrightarrow d_i \in (0, \epsilon]$.

We now give the overall algorithm which is based on a separation into boxes that are d - δ -big in all directions and boxes that are d - δ -small in at least one direction. We explain the details in the proof only.

Algorithm E

1. Set $\delta := \epsilon/\lceil 37(7 + \epsilon) \rceil$, let

$$L_1 := \{R_i \mid R_i \text{ is } d\text{-}\delta\text{-big for each } d \in \{x, y, z\}\}$$

and $L_2 := L \setminus L_1$.

2. For each $L_3 \subseteq L_1$ such that $|L_3| \leq \lfloor 1/\delta^3 \rfloor$ uses an exact algorithm to test the packability of L_3 . Store a feasible L_3 of maximum total profit.
3. Use an FPTAS for classical KP from [21, 27] to select $L_4 \subseteq L_2$ such that $V(L_4) \leq 1$, and $P(L_4) \geq (1 - \delta)OPT$ holds.
4. Use the construction described below to select $L_5 \subseteq L_4$ which can be packed into at most 6 bins at a small loss of profit.
5. Out of the at most 7 sets generated in Step 2, and Step 4 returns one with the maximum profit.

Theorem 13. *Algorithm E is a $(7 + \epsilon)$ -approximation algorithm for OKP-3.*

Proof. Note that $\lfloor 1/\delta^3 \rfloor$ is an upper bound for the number of boxes from L_1 in a feasible solution since δ^3 is a lower bound for the volume of each $R_i \in L_1$. Step 2 can be carried out in time polynomial in δ and thus polynomial in $1/\epsilon$ using an exact optimization algorithm as in [18]. We show that in Step 4 at most 6 sets

are generated, resulting in at most 7 bins in total. Partition L_4 into 3 subsets X' , Y' and Z' such that in each of these all boxes R_i are d - δ -small for the corresponding direction; note that $V(X') + V(Y') + V(Z') \leq 1$ holds. We apply the construction from Theorem 12 in each of the three directions. Study the following cases, where $V(X') \geq V(Y') \geq V(Z')$ holds w.l.o.g.

Case 1. $V(X') \in (3/4, 1]$. The boxes in X' can be packed into at most 4 bins. We have $V(Y') + V(Z') < 1/4$. This means $V(Y') < 1/4$ and $V(Z') < 1/4$ hold. Consequently Y' and Z' can be packed into at most one bin each, resulting in at most 7 bins.

Case 2. $V(X') \in (1/2, 3/4]$. The boxes in X' can be packed into at most three bins. Furthermore, $V(Y') + V(Z') < 1/2$, which means that $V(Y') < 1/2$ holds. Consequently the boxes in Y' can be packed into at most 2 bins. Furthermore, $V(Z') < 1/4$ holds and finally the boxes in Z' can be packed into at most 1 bin; this generates at most 7 bins in total.

Case 3. We have $V(X') \in [0, 1/2]$. The boxes in X' can be packed into at most two additional bins. Furthermore, $V(Y') \leq 1/2$ and $V(Z') \leq 1/2$ hold. This means that the boxes in Y' and Z' can be packed into at most two bins each. In total at most 7 bins are generated.

To prove the approximation ratio we fix an optimal solution S , and let P_1^* be the profit of boxes in $S \cap L_1$ and P_2^* be the profit of boxes in $S \cap L_2$. Consequently $P_1^* + P_2^* = OPT$ holds. Let P_1 be the profit of the set that is stored in Step 2 and let P_2 be the profit of the set that is selected in Step 3. By construction we have $P_1 \geq P_1^*$ and $P_2 \geq (1 - \delta)P_2^*$. Furthermore, by three-fold applications of the construction from Theorem 12 the loss of profit in P_2 is bounded by $36\delta P_2$; let P_2' denote the remaining list. The profit of the set returned in Step 5 is at least

$$\begin{aligned} \frac{1}{7}(P_1 + P_2') &\geq \frac{1}{7}(P_1^* + (1 - \delta)(1 - 36\delta)P_2^*) \\ &\geq \frac{1}{7}(P_1^* + P_2^*)(1 - \delta)(1 - 36\delta) \\ &= \frac{1}{7}OPT(1 - \delta)(1 - 36\delta) \geq \frac{1}{(7 + \epsilon)}OPT, \end{aligned}$$

which proves the claimed approximation ratio. \square

Theorem 14. *The bound of $(7 + \epsilon)$ for the approximation ratio of Algorithm E is asymptotically tight in the sense that it cannot be improved for ϵ arbitrary small.*

Proof. Let $\epsilon \in (0, 1/4]$ and δ be defined as in Step 1 of Algorithm E. Note that $\delta \leq 1/259$. Set $\gamma := \max\{(12i)^{-1} \mid i \in \mathbb{N}, (12i)^{-1} \leq \delta\}$. Let $\alpha \in \mathbb{R}_+$ such that $(3/4 - \alpha)(1/3 + \alpha) > 1/4$, $(1/3 - 2\alpha)(3/4 + 6\alpha) >$

$1/4$ and $2\alpha \leq \gamma$ hold. It is easy to see that such an α exists. Note that $\{1/(2\gamma), 1/(3\gamma), 1/(4\gamma), 1/(6\gamma)\} \subseteq \mathbb{N}$. We use the boxes,

$$\begin{aligned} A_i &:= \left(\frac{1}{2}, \frac{1}{2} + \alpha, \gamma\right), \quad i \in \left\{1, \dots, \frac{1}{\gamma}\right\}, \\ A &:= \{A_1, \dots, A_{\frac{1}{\gamma}}\}, \\ B_i &:= \left(\frac{1}{2}, \frac{1}{2} + \alpha, \gamma\right), \quad i \in \left\{1, \dots, \frac{1}{4\gamma} - 1\right\}, \\ B &:= \{B_1, \dots, B_{\frac{1}{4\gamma} - 1}\}, \\ C_i &:= \left(\gamma, \frac{1}{2} + \alpha, \frac{1}{2}\right), \quad i \in \left\{1, \dots, \frac{1}{2\gamma}\right\}, \\ C &:= \{C_1, \dots, C_{\frac{1}{2\gamma}}\}, \\ D_i &:= \left(\gamma, 1, \frac{1}{4} + \alpha\right), \quad i \in \left\{1, \dots, \frac{1}{2\gamma}\right\}, \\ D &:= \{D_1, \dots, D_{\frac{1}{2\gamma}}\}, \\ E_i &:= \left(\gamma, \frac{1}{3} + \alpha, \frac{3}{4} - \alpha\right), \quad i \in \left\{1, \dots, \frac{1}{3\gamma}\right\}, \\ E &:= \{E_1, \dots, E_{\frac{1}{3\gamma}}\}, \\ F_i &:= \left(\frac{1}{3} + \alpha, \gamma, \frac{3}{4} - \alpha\right), \quad i \in \left\{1, \dots, \frac{1}{6\gamma} - 1\right\}, \\ F &:= \{F_1, \dots, F_{\frac{1}{6\gamma} - 1}\}, \\ G_i &:= \left(\frac{1}{3} + \alpha, \gamma, \frac{3}{4} - \alpha\right), \quad i \in \left\{1, \dots, \frac{1}{2\gamma} - 1\right\}, \\ G &:= \{G_1, \dots, G_{\frac{1}{2\gamma} - 1}\}, \\ H_i &:= \left(\frac{1}{3} - 2\alpha, \gamma, \frac{3}{4} + 6\alpha\right), \quad i \in \left\{1, \dots, \frac{1}{2\gamma} - 1\right\}, \\ H &:= \{H_1, \dots, H_{\frac{1}{2\gamma} - 1}\}, \\ R_1 &:= (\delta + \alpha, \delta + \alpha, \delta + \alpha), \end{aligned}$$

to define the list L . In the following we will show that L admits a feasible packing into the unit cube. Note that each of the sets (without the box R_1) defined above can be arranged to a rectangular block by stacking them on top of each other in the direction d in which they are d - δ -small, see Fig.6. Furthermore, we note that

$$\begin{aligned} |A| &= \frac{1}{\gamma}, br_z(A_i) > \frac{1}{4} \quad \text{for } A_i \in A, \sum_{A_i \in A} z(A_i) = 1, \\ |B| &= \frac{1}{4\gamma} - 1, br_z(B_i) > \frac{1}{4} \quad \text{for } B_i \in B, \\ &\sum_{B_i \in B} z(B_i) = \frac{1}{4} - \gamma, \\ |C| &= \frac{1}{2\gamma}, br_x(C_i) > \frac{1}{4} \quad \text{for } C_i \in C, \sum_{C_i \in C} x(C_i) = \frac{1}{2}, \end{aligned}$$

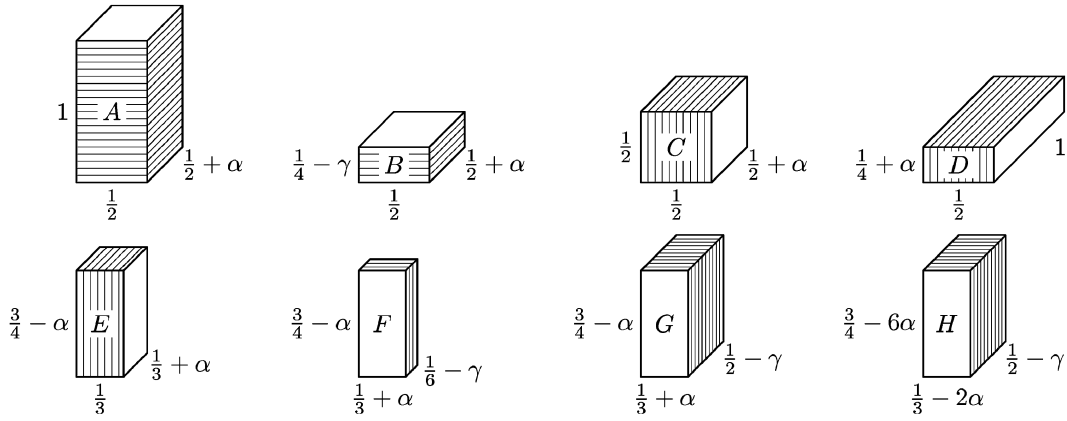


Fig.6. Boxes arranged in blocks.

$$\begin{aligned}
 |D| &= \frac{1}{2\gamma}, br_x(D_i) > \frac{1}{4} \quad \text{for } D_i \in D, \sum_{D_i \in D} x(D_i) = \frac{1}{2}, \\
 |E| &= \frac{1}{3\gamma}, br_x(E_i) > \frac{1}{4} \quad \text{for } E_i \in E, \sum_{E_i \in E} x(E_i) = \frac{1}{3}, \\
 |F| &= \frac{1}{6\gamma} - 1, br_y(F_i) > \frac{1}{4} \quad \text{for } F_i \in F, \\
 &\sum_{F_i \in F} y(F_i) = \frac{1}{6} - \gamma, \\
 |G| &= \frac{1}{2\gamma} - 1, br_y(G_i) > \frac{1}{4} \quad \text{for } G_i \in G, \\
 &\sum_{G_i \in G} y(G_i) = \frac{1}{2} - \gamma, \\
 |H| &= \frac{1}{2\gamma} - 1, br_y(H_i) > \frac{1}{4} \quad \text{for } H_i \in H, \\
 &\sum_{H_i \in H} y(H_i) = \frac{1}{2} - \gamma
 \end{aligned}$$

hold, where $br_d(R)$ denotes the base rectangle in direction d and $d(R)$ denotes component d for a box R , and $d \in \{x, y, z\}$. As indicated above, these sets together with the box R_1 permit a feasible packing since

- F, E, C and F, E, B can be placed next to one another in direction y ,
- C, B, D can be placed next to one another in direction z ,
- A, H and A, G and B, G and C, G can be placed next to one another in direction x ,

which can be verified by elementary calculation, see Fig.7.

In the following we discuss the execution of Algorithm E. The profits necessary for our construction will be described together with the presentation. First we

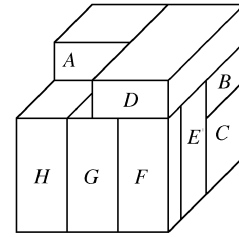


Fig.7. Feasible packing of the boxes into the unit cube.

require $P(L) = 1$; this also equals OPT since L permits a feasible packing. In Step 1 we obtain $L_1 = \{R_1\}$ and $L_2 = L \setminus \{R_1\}$. Step 2 stores $\{R_1\}$ as a feasible candidate for selection later; here we require $P(R_1) = 1/(7 + \epsilon)$. In Step 3 we assume that $L_4 = L_2$ which is possible since L_2 is an optimal solution for the generated knapsack instance; note that here

$$P(L_4) = 1 - 1/(7 + \epsilon) = (6 + \epsilon)/(7 + \epsilon).$$

Now we discuss how L_4 is packed in Step 4. First L_4 is partitioned into

$$X' = C \cup D \cup E, \quad Y' = F \cup G \cup H, \quad Z' = A \cup B.$$

Then each of these three sets is packed into a strip in the corresponding direction. Note that

$$\begin{aligned}
 br_x(R) &> \frac{1}{4} \quad \text{for } R \in X', \\
 br_y(R) &> \frac{1}{4} \quad \text{for } R \in Y', \\
 br_z(R) &> \frac{1}{4} \quad \text{for } R \in Z'
 \end{aligned}$$

hold. Hence Algorithm E packs all boxes on top of one another in the corresponding direction; we assume that the boxes in each of the sets are packed in “lexicographical” order. Note that the first layer is separated. We denote the heights of the generated strips for X' , Y'

and Z' by $h_{X'}$, $h_{Y'}$ and $h_{Z'}$, respectively, and obtain

$$h_{X'} = \frac{4}{3} - \gamma, \quad h_{Y'} = \frac{7}{6} - 4\gamma, \quad h_{Z'} = \frac{5}{4} - 2\gamma.$$

Since $\gamma \leq \delta \leq 1/259$, all of these strips have height more than 1 and will later result in 2 bins. Let X'_1 denote the set of boxes that are in or intersect the region $[0, 1] \times [0, 1] \times [0, 1]$ in the strip and set $X'_2 := X' \setminus X'_1$. We define Y'_1 , Y'_2 and Z'_1 , Z'_2 in a similar way. Now we require that

$$\begin{aligned} P(X'_1) &= P(X'_2) = P(Y'_1) = P(Y'_2) \\ &= P(Z'_1) = P(Z'_2) = \frac{1}{6} \frac{6 + \epsilon}{7 + \epsilon}. \end{aligned}$$

The profit is evenly partitioned among the items in the sets. Algorithm E uses the shifting technique from Lemma 11 to generate feasible packings for X' , Y' and Z' , respectively, into 2 bins. Hence three gaps of height δ are generated into each of the strips causing a small loss of profit; the separated layer and the boxes intersecting the cutting square are merged into each strip. It is easy to see that the shifting technique acts only on X'_1 , Y'_1 , and Z'_1 , respectively, since the profit on each items here is lower than in the corresponding second sets. Thus no item is swapped between sets by the shifting technique. We obtain 6 bins, each containing subsets of X'_1 , X'_2 , Y'_1 , Y'_2 , Z'_1 , and Z'_2 .

Since each of the sets holds a profit of $(6 + \epsilon)/[6(7 + \epsilon)]$, each bin holds a profit less than or equal to $(1 + \epsilon/6)/(7 + \epsilon)$ which finishes the proof. \square

7 Rotational Case

Finally, we discuss the application of Algorithm E in two different rotational scenarios. In both scenarios rotations of the boxes are only permitted by 90° around certain axes. In the first case, which we denote by *z-oriented* OKP-3, rotations are only permitted around the z -axis. This setting is motivated by packing fragile goods and has been considered in the strip packing variant in [29–31]. In the second case, which we denote by *rotational* OKP-3, rotations are permitted around all three axes. Note that for both scenarios $Q = (1, 1, 1)$ does not hold without loss of generality, but is an explicit assumption. Surprisingly, a better approximation ratio can be obtained easily although implicitly, the search space is dramatically enlarged.

We show how Algorithm E can be revised to yield a better approximation ratio for both scenarios. Step 2 is modified in such a way that the exact packing algorithm takes rotations into account, and as before, here

at most 1 bin is generated. The most important part is the modification of Step 4. We separate the description of the two scenarios and start with the *z-oriented* setting. Let Z be the set of boxes $R_i \in L_4$ that are z - δ -small and $X = L_4 \setminus Z$. We introduce a preprocessing step in which each $R_i \in X$ is rotated in such a way that the side length x_i is minimal. Consequently, $x_i \leq \delta$ holds for each $R_i \in X$. Hence, the generation of a strip for direction y can be removed; we build only two strips in directions x and z to which the shifting technique from Theorem 12 is applied. This again causes an additional loss of profit which is bounded, however. Let $i \in \{1, \dots, 4\}$ such that $(i - 1)/4 \leq V(Z) \leq i/4$. Then by Theorem 12, exactly i bins are needed to pack the strip in direction z . Furthermore, we get

$$V(X) \leq V(L_4) - V(Z) \leq 1 - \frac{i - 1}{4} = \frac{5 - i}{4},$$

which yields a number of $5 - i$ bins for the corresponding strip. In total a number of $1 + i + (5 - i) = 6$ bins are needed to pack all items. We call the resulting approach Algorithm F and obtain the following result.

Theorem 15. *Algorithm F is a $(6 + \epsilon)$ -approximation algorithm for *z-oriented* OKP-3.*

For the *rotational* OKP-3 we have to continue the approach above. Since rotations around all axes are permitted, all items that are d - δ -small for any direction $d \in \{x, y, z\}$ can be packed into one strip in direction z by orienting them such that they are z - δ -small. Since $V(L_4) \leq 1$, application of the construction of Theorem 12 yields at most 4 additional bins. We call the resulting approach Algorithm G and obtain the following result.

Theorem 16. *Algorithm G is a $(5 + \epsilon)$ -approximation algorithm for *rotational* OKP-3.*

8 Conclusion

We contributed approximation algorithms for an NP-hard combinatorial optimization problem, where the runtimes of the simpler algorithms are practical. It is of interest whether here an algorithm with ratio $6 + \epsilon$ or less exists. We are interested in a reduction of the running time, especially for Algorithm E. In [32] it was proved that it is NP-complete to decide whether a set of squares can be packed into the unit square. However, it is an open problem whether checking the packability of cubes into the unit cube is NP-complete.

Acknowledgements The authors thank the anonymous referees for valuable comments. Florian Diedrich thanks Denis Naddef for hospitality during the preparation of this article.

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